

Interference Minimization in Asymmetric Sensor Networks^{*}

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Abstract. A fundamental problem in wireless sensor networks is to connect a given set of sensors while minimizing the *receiver interference*. This is modeled as follows: each sensor node corresponds to a point in \mathbb{R}^d and each *transmission range* corresponds to a ball. The receiver interference of a sensor node is defined as the number of transmission ranges it lies in. Our goal is to choose transmission radii that minimize the maximum interference while maintaining a strongly connected asymmetric communication graph.

For the two-dimensional case, we show that it is NP-complete to decide whether one can achieve a receiver interference of at most 5. In the one-dimensional case, we prove that there are optimal solutions with non-trivial structural properties. These properties can be exploited to obtain an exact algorithm that runs in quasi-polynomial time. This generalizes a result by Tan et al. to the asymmetric case.

1 Introduction

Wireless sensor networks constitute a popular paradigm in mobile networks: several small independent devices are distributed in a certain region, and each device has limited computational resources. The devices can communicate through a wireless network. Since battery life is limited, it is imperative that the overhead for the communication be kept as small as possible. One major concern when trying to achieve this goal is to control the *interference* caused by competing senders. This enables us to reduce the range of the senders, thus increasing battery life. At the same time, we need to ensure that the resulting communication graph remains connected.

There are many different ways to formalize the problem of interference minimization. Usually, the devices are modeled as points in d -dimensional space, and the transmission ranges are modeled as d -dimensional balls. Each point can choose the radius of its transmission range, and different choices of transmission

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ranges lead to different reachability structures. There are two ways to interpret the resulting communication graph. In the *symmetric* case, the communication graph is undirected, and it contains an edge between two points p and q if and only if both p and q lie in the transmission range of the other point. For a valid assignment of transmission ranges, we require that the communication graph is connected. In the *asymmetric* case, the communication graph is directed, and there is an edge from p to q if and only if p lies in the transmission range of q . We require that the communication graph is strongly connected, or, in a slightly different model, that there is one point that is reachable from every other point through a directed path.

In both the symmetric and the asymmetric case, the (*receiver-centric*) *interference* of a point is defined as the number of transmission ranges that it lies in [5]. The goal is to find a valid assignment of transmission ranges that makes the maximum interference as small as possible. We refer to the resulting interference as *minimum interference*. The minimum interference under the two models for the asymmetric case differs by at most one: if there is a point reachable from every other, we can increase its transmission range to include all other points. As a result, the communication graph becomes strongly connected, while the minimum interference increases by at most one.

Let n be the number of points. In the symmetric case, one can always achieve interference $O(\sqrt{n})$, and this is sometimes necessary [2, 5]. In the one-dimensional case, there is an efficient approximation algorithm with approximation factor $O(n^{1/4})$ [5]. Furthermore, Tan et al. [6] prove the existence of optimal solutions with interesting structural properties in one dimension. This can be used to obtain a nontrivial exact algorithm for this case. In the asymmetric case, the interference is significantly smaller: one can always achieve interference $O(\log n)$, which is sometimes optimal (e.g., [3]).

Our results. We consider interference minimization in asymmetric wireless sensor networks in one and two dimensions. We show that for two dimensions, it is NP-complete to find a valid assignment that minimizes the maximum interference. In one dimension we consider our second model requiring one point that is reachable from every other point through a directed path. Generalizing the result by Tan et al. [6], we show that there is an optimal solution that exhibits a certain binary tree structure. By means of dynamic programming, this structure can be leveraged for a nontrivial exact algorithm. Unlike the symmetric case, this algorithm always runs in quasi-polynomial time $2^{O(\log^2 n)}$, making it unlikely that the one-dimensional problem is NP-hard. Nonetheless, a polynomial time algorithm remains elusive.

2 Preliminaries and Notation

We now formalize our interference model for the planar case. Let $P \subset \mathbb{R}^2$ be a planar n -point set. A *receiver assignment* $N : P \rightarrow P$ is a function that assigns to each point in P the furthest point that receives data from P . The resulting (asymmetric) *communication graph* $G_P(N)$ is the directed graph with

vertex set P and edge set $E_P(N) = \{(p, q) \mid \|p - q\| \leq \|p - N(p)\|\}$, i.e., from each point $p \in P$ there are edges to all points that are at least as close as the assigned receiver $N(p)$. The receiver assignment N is *valid* if $G_P(N)$ is strongly connected.

For $p \in \mathbb{R}^2$ and $r > 0$, let $B(p, r)$ denote the closed disk with center p and radius r . We define $B_P(N) = \{B(p, d(p, N(p))) \mid p \in P\}$ as the set that contains for each $p \in P$ a disk with center p and $N(p)$ on the boundary. The disks in $B_P(N)$ are called the *transmission ranges* for N . The *interference* of N , $I(N)$, is the maximum number of transmission ranges that cover a point in P , i.e., $I(N) = \max_{p \in P} |\{p \in B \mid B \in B_P(N)\}|$. In the *interference minimization problem*, we are looking for a valid receiver assignment with minimum interference.

3 NP-completeness in Two Dimensions

We show that the following problem is NP-complete: given a planar point set P , does there exist a valid receiver assignment N for P with $I(N) \leq 5$? It follows that the minimum interference for planar point sets is NP-hard to approximate within a factor of $6/5$.

The problem is clearly in NP. To show that interference minimization is NP-hard, we reduce from the problem of deciding whether a grid graph of maximum degree 3 contains a Hamiltonian path: a *grid graph* G is a graph whose vertex set $V \subset \mathbb{Z} \times \mathbb{Z}$ is a finite subset of the integer grid. Two vertices $u, v \in V$ are adjacent in G if and only if $\|u - v\|_1 = 1$, i.e., if u and v are neighbors in the integer grid. A *Hamiltonian path* in G is a path that visits every vertex in V exactly once. Papadimitriou and Vazirani showed that it is NP-complete to decide whether a grid graph G of maximum degree 3 contains a Hamiltonian cycle [4]. Note that we may assume that G is connected; otherwise there can be no Hamiltonian path.

Our reduction proceeds by replacing each vertex v of the given grid graph G by a *vertex gadget* P_v ; see Fig. 1. The vertex gadget consists of 13 points, and it has five parts: (a) the *main point* M with the same coordinates as v ; (b) three

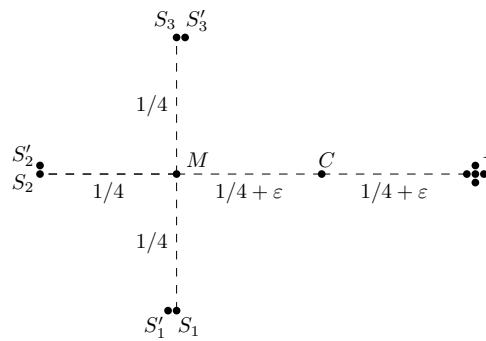


Fig. 1. The vertex gadget.

satellite stations with two points each: $S_1, S'_1, S_2, S'_2, S_3, S'_3$. The coordinates of the S_i are chosen from $\{v \pm (0, 1/4), v \pm (1/4, 0)\}$ so that there is a satellite station for each edge in G that is incident to v . If v has degree two, the third satellite station can be placed in any of the two remaining directions. The S'_i lie at the corresponding clockwise positions from $\{v \pm (\varepsilon, 1/4), v \pm (1/4, -\varepsilon)\}$, for a sufficiently small $\varepsilon > 0$; (c) the *connector* C , a point that lies roughly at the remaining position from $\{v \pm (0, 1/4), v \pm (1/4, 0)\}$ that is not occupied by a satellite station, but an ε -unit further away from M . For example, if $v + (0, 1/4)$ has no satellite station, then C lies at $v + (0, 1/4 + \varepsilon)$; and (d) the *inhibitor*, consisting of five points I_c, I_1, \dots, I_4 . The point I_c is the center of the inhibitor and I_1 is the point closest to C . The position of I_c is $M + 2(C - M) + \varepsilon(C - M)/\|C - M\|$, that is, the distance between I_c and C is an ε -unit larger than the distance between C and M : $\|M - C\| + \varepsilon = \|C - I_c\|$. The points I_1, \dots, I_4 are placed at the positions $\{I_c \pm (0, \varepsilon), I_c \pm (\varepsilon, 0)\}$, with I_1 closest to C .

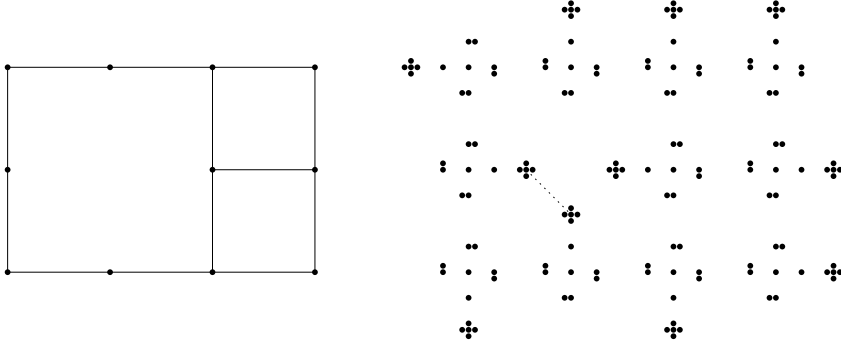


Fig. 2. An example reduction.

Given a grid graph G , the reduction can be carried out in polynomial time: just replace each vertex v of G by the corresponding gadget P_v ; see Fig. 2 for an example. Let $P = \bigcup_{v \in G} P_v$ be the resulting point set. Two satellite stations in P that correspond to the same edge of G are called *partners*. First, we investigate the interference in any valid receiver assignment for P .

Lemma 3.1. *Let N be a valid receiver assignment for P . Then in each vertex gadget, the points I_c and M have interference as least 5, and the points S_1, S_2 , and S_3 have interference at least 3.*

Proof. For each point $p \in P$, the transmission range $B(p, d(p, N(p)))$ must contain at least the nearest neighbor of p . Furthermore, in each satellite station and in each inhibitor, at least one point must have an assigned receiver outside of the satellite station or inhibitor; otherwise, the communication graph $G_P(N)$ would not be strongly connected. This forces interference of 5 at M and at I_c : each satellite station and C must have an edge to M , and I_1, \dots, I_4 all must have an

edge to I_c . Similarly, for $i = 1, \dots, 3$, the main point M and the satellite S'_i must have an edge to S_i ; see Fig. 3. \square

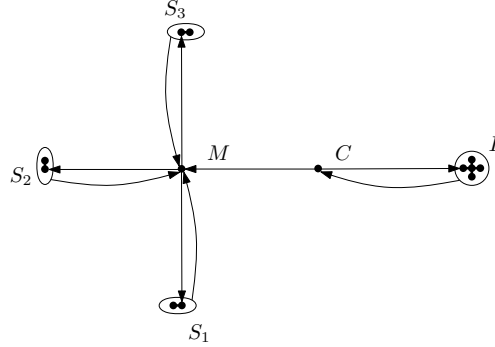


Fig. 3. The nearest neighbors in a vertex gadget.

Let N be a valid receiver assignment, and let P_v be a vertex gadget in P . An *outgoing* edge for P_v is an edge in $G_P(N)$ that originates in P_v and ends in a different vertex gadget. An *incoming* edge for P_v is an edge that originates in a different gadget and ends in P_v . A *connecting* edge for P_v is either an outgoing or an incoming edge for P_v . If $I(N) \leq 5$ holds, then Lemma 3.1 implies that a connecting edge can be incident only to satellite stations. The proof of the following lemma is given in Appendix A.

Lemma 3.2. *Let N be a valid receiver assignment for P with $I(N) \leq 5$. Let P_v be a vertex gadget of P and e an outgoing edge from P_v to another vertex gadget P_w . Then e goes from a satellite station of P_v to its partner satellite station in P_w . Furthermore, in each satellite station of P_v , at most one point is incident to outgoing edges.*

Next, we show that the edges between the vertex gadgets are quite restricted.

Lemma 3.3. *Let N be a valid receiver assignment for P with $I(N) \leq 5$. For every vertex gadget P_v in P , at most two satellite stations in P_v are incident to connecting edges in $G_P(N)$.*

Proof. By Lemma 3.2 connecting edges are between satellite stations and by Lemma 3.1, the satellite points S_i in P_v have interference at least 3.

First, assume that all three satellite stations in P_v have outgoing edges. This would increase the interference at all three S_i to 5. Then, P_v could not have any incoming edge from another vertex gadget, because this would increase the interference for at least one S_i (note that due to the placement of the S'_i , every incoming edge causes interference at an S_i). If P_v had no incoming edge, $G_P(N)$

would not be strongly connected. It follows that P_v has at most two satellite stations with outgoing edges.

Next, assume that two satellite stations in P_v have outgoing edges. Then, the third satellite station of P_v cannot have an incoming edge, as the two outgoing edges already increase the interference at the third satellite station to 5.

Hence, we know that every vertex gadget P_v either (i) has connecting edges with all three partner gadgets, exactly one of which is outgoing, or (ii) is connected to at most two other vertex gadgets. Take a vertex gadget P_v of type (i) with partners P_{u_1}, P_{u_2}, P_w . Suppose that P_v has incoming edges from P_{u_1} and P_{u_2} and that the outgoing edge goes to P_w . Follow the outgoing edge to P_w . If P_w is of type (i), follow the outgoing edge from P_w ; if P_w is of type (ii) and has an outgoing edge to a vertex gadget we have not seen yet, follow this edge. Continue this process until P_v is reached again or until the next vertex gadget has been visited already. This gives all vertex gadgets that are reachable from P_v on a directed path. However, in each step there is only one choice for the next vertex gadget. Thus, the process cannot discover P_{u_1} and P_{u_2} , since both of them would lead to P_v in the next step, causing the process to stop. It follows that at least one of P_{u_1} or P_{u_2} is not reachable from P_v , although $G_P(N)$ should be strongly connected. Therefore, all vertex gadgets in $G_P(N)$ must be of type (ii), as claimed in the lemma. \square

We can now prove the main theorem of this section.

Theorem 3.4. *Given a point set $P \subset \mathbb{R}^2$, it is NP-complete to decide whether there exists a valid receiver assignment N for P with $I(N) \leq 5$.*

Proof. Using the receiver assignment N as certificate, the problem is easily seen to be in NP. To show NP-hardness, we use the polynomial time reduction from the Hamiltonian path problem in grid graphs: given a grid graph G of maximum degree 3, we construct a planar point set P as above. It remains to verify that G has a Hamiltonian path if and only if P has a valid receiver assignment N with $I(N) \leq 5$.

Given a Hamilton path H in G , we construct a valid receiver assignment N for P as follows: in each vertex gadget, we set $N(M) = C$, $N(C) = M$, and $N(I_1) = C$. For $i = 1, \dots, 3$ we set $N(S'_i) = S_i$ and $N(I_{i+1}) = I_c$. Finally, we set $N(I_c) = I_1$. This essentially creates the edges from Fig. 3, plus the edge from M to C . Next, we encode H into N : for each S_i on an edge of H , we set $N(S_i)$ to the corresponding S_i in the partner station. For the remaining S_i , we set $N(S_i) = M$. Since H is Hamiltonian, $G_P(N)$ is strongly connected (note that each vertex gadget induces a strongly connected subgraph). It can now be verified that M and I_c have interference 5; I_2, I_3, I_4 have interference 2; and I_1 has interference 3. The point C has interference between 2 and 4, depending on whether S_1 and S_3 are on edges of H . The satellites S_i and S'_i have interference at most 5 and 4, respectively.

Now consider a valid receiver assignment N for P with $I(N) \leq 5$. Let F be the set of edges in G that correspond to pairs of vertex gadgets with a connecting edge in $G_P(N)$. Let H be the subgraph that F induces in G . By Lemma 3.3,

H has maximum degree 2. Furthermore, since $G_P(N)$ is strongly connected, the graph H is connected and meets all vertices of G . Thus, H is a Hamiltonian path (or cycle) for G , as desired. \square

Remark. A similar result to Theorem 3.4 also holds for symmetric communication graphs networks [1].

4 The One-Dimensional Case

For the one-dimensional case we minimize receiver interference under the second model discussed in the introduction: given $P \subset \mathbb{R}$ and a receiver assignment $N : P \rightarrow P$, the graph $G_P(N)$ now has a directed edge from each point $p \in P$ to its assigned receiver $N(p)$, and no other edges. N is *valid* if $G_P(N)$ is acyclic and if there is a sink $r \in P$ that is reachable from every point in P . The sink has no outgoing edge. The interference of N , $I(N)$, is defined as before.

4.1 Properties of Optimal Solutions

We now explore the structure of optimal receiver assignments. Let $P \subset \mathbb{R}$ and N be a valid receiver assignment for P with sink r . We can interpret $G_P(N)$ as a directed tree, so we call r the *root* of $G_P(N)$. For a directed edge pq in $G_P(N)$, we say that p is a *child* of q and q is the *parent* of p . We write $p \rightsquigarrow_N q$ if there is a directed path from p to q in $G_P(N)$. If $p \rightsquigarrow_N q$, then q is an *ancestor* of p and p a *descendant* of q . Note that p is both an ancestor and a descendant of p . Two points $p, q \in P$ are *unrelated* if p is neither an ancestor nor a descendant of q . For two points p, q , we define $((p, q)) = (\min\{p, q\}, \max\{p, q\})$ as the open interval bounded by p and q , and $[[p, q]] = [\min\{p, q\}, \max\{p, q\}]$ as the closure of $((p, q))$. An edge pq of $G_P(N)$ is a *cross edge* if the interval $((p, q))$ contains at least one point that is not a descendant of p .

Our main structural result is that there is always an optimal receiver assignment for P without cross edges. A similar property was observed by Tan et al. for the symmetric case [6].

Lemma 4.1. *Let N^* be a valid receiver assignment for P with minimum interference. There is a valid assignment \tilde{N} for P with $I(\tilde{N}) = I(N^*)$ such that $G_P(\tilde{N})$ has no cross edges.*

Proof. Pick a valid assignment \tilde{N} with minimum interference that minimizes the total length of the cross edges

$$C(\tilde{N}) := \sum_{pq \in C} \|p - q\|,$$

where C are the cross-edges of $G_P(\tilde{N})$. If $C(\tilde{N}) = 0$, we are done. Thus, suppose $C(\tilde{N}) > 0$. Pick a cross edge pq such that the hop-distance (i.e., the number of edges) from p to the root is maximum among all cross edges. Let p_l be the leftmost and p_r the rightmost descendant of p . We refer to Appendix A for a proof of the following lemma.

Proposition 4.2 *The interval $[p_l, p_r]$ contains only descendants of p .*

Let R be the points in $((p, q))$ that are not descendants of p . Each point in R is either unrelated to p , or it is an ancestor of p . Let $z \in R$ be the point in R that is closest to p (i.e., z either lies directly to the left of p_l or directly to the right of p_r). We now describe how to construct a new valid assignment \hat{N} , from which we will eventually derive a contradiction to the choice of \tilde{N} . The construction is as follows: replace the edge pq by pz . Furthermore, if (i) $q \rightsquigarrow_{\tilde{N}} z$; (ii) the last edge $z'z$ on the path from q to z crosses the interval $[p_l, p_r]$; and (iii) $z'z$ is not a cross-edge, we also change the edge $z'z$ to the edge that connects z' to the closer of p_l or p_r . We give the proof of the following proposition in Appendix A.

Proposition 4.3 *\hat{N} is a valid assignment.*

Proposition 4.4 *We have $I(N^*) = I(\hat{N})$.*

Proof. Since the new edges are shorter than the edges they replace, each transmission range for \hat{N} is contained in the corresponding transmission range for \tilde{N} . The interference cannot decrease since N^* is optimal. \square

Proposition 4.5 *We have $C(\hat{N}) < C(\tilde{N})$.*

Proof. First, we claim that \hat{N} contains no new cross edges, except possibly pz : suppose ab is a cross edge of $G_P(\hat{N})$, but not of $G_P(\tilde{N})$. This means that $((a, b))$ contains a point x with $x \rightsquigarrow_{\tilde{N}} a$, but $x \not\rightsquigarrow_{\hat{N}} a$. Then x must be a descendant of p in $G_P(\tilde{N})$ and in $G_P(\hat{N})$, because as we saw in the proof of Claim 4.3, for any $y \in P \setminus [p_l, p_r]$, we have that if $y \rightsquigarrow_{\tilde{N}} a$, then $y \rightsquigarrow_{\hat{N}} a$.

Hence, $((a, b))$ and $[p_l, p_r]$ intersect. Since ab is a cross edge, the choice of pq now implies that $[p_l, p_r] \subseteq ((a, b))$. Thus, z lies in $[[a, b]]$, because z is a direct neighbor of p_l or p_r . We claim that $b = z$. Indeed, otherwise we would have $z \rightsquigarrow_{\tilde{N}} a$ (since ab is not a cross edge in $G_P(\tilde{N})$), and thus also $z \rightsquigarrow_{\hat{N}} a$. However, we already observed $x \rightsquigarrow_{\hat{N}} p$, so we would have $x \rightsquigarrow_{\hat{N}} a$ (recall that we introduce the edge pz in \hat{N}). This contradicts our choice of x .

Now it follows that $ab = az$ is the last edge on the path from p to z , because if a were not an ancestor of p , then ab would already be a cross-edge in $G_P(\tilde{N})$. Hence, (i) a is an ancestor of q ; (ii) az crosses the interval $[p_l, p_r]$; and (iii) az is not a cross edge in \tilde{N} . These are the conditions for the edge $z'z$ that we remove from \tilde{N} . The new edge e from a to p_l or p_r cannot be a cross edge, because ab is not a cross edge in $G_P(\hat{N})$ and e does not cover any descendants of p .

Hence, $G_P(\hat{N})$ contain no new cross-edges, except possibly pz which replaces the cross edge pq . By construction, $\|p - z\| < \|p - q\|$, so $C(\hat{N}) < C(\tilde{N})$. \square

Propositions 4.3–4.5 yield a contradiction to the choice of \tilde{N} . It follows that we must have $C(\tilde{N}) = 0$, as desired. \square

Let $P \subset \mathbb{R}$. We say that a valid assignment N for P has the *BST-property* if the the following holds for any vertex p of $G_P(N)$: (i) p has at most one child q

with $p < q$ and at most one child q with $p > q$; and (ii) let p_l be the leftmost and p_r the rightmost descendant of p . Then $[p_l, p_r]$ contains only descendants of p . In other words: $G_P(N)$ constitutes a binary search tree for the (coordinates of the) points in P . A valid assignment without cross edges has the BST-property. The following is therefore an immediate consequence of Lemma 4.1.

Theorem 4.6. *Every $P \subset \mathbb{R}$ has an optimal valid assignment with the BST-property.* \square

4.2 A Quasi-Polynomial Algorithm

We now show how to use Theorem 4.6 for a quasi-polynomial time algorithm to minimize the interference. The algorithm uses dynamic programming. A subproblem π for the dynamic program consists of four parts: (i) an interval $P_\pi \subseteq P$ of *consecutive* points in P ; (ii) a root $r_\pi \in P_\pi$; (iii) a set I_π of *incoming interference*; and (iv) a set O_π of *outgoing interference*.

The objective of π is to find an optimal valid assignment N for P_π subject to (i) the root of $G_N(P_\pi)$ is r_π ; (ii) the set O_π contains all transmission ranges of $B_{P_\pi}(N)$ that cover points in $P \setminus P_\pi$ plus potentially a transmission range with center r_π ; (iii) the set I_π contains transmission ranges that cover points in P_π and have their center in $P \setminus P_\pi$. The interference of N is defined as the maximum number of transmission ranges in $B_{P_\pi}(N) \cup I_\pi \cup O_\pi$ that cover any given point of P_π . The transmission ranges in $O_\pi \cup I_\pi$ are given as pairs $(p, q) \in P^2$, where p is the center and q a point on the boundary of the range.

Each range in $O_\pi \cup I_\pi$ covers a boundary point of P_π . Since it is known that there is always an assignment with interference $O(\log n)$ (see [5] and Observation B.1), no point of P lies in more than $O(\log n)$ ranges of $B_P(N^*)$. Thus, we can assume that $|I_\pi \cup O_\pi| = O(\log n)$, and the total number of subproblems is $n^{O(\log n)}$.

A subproblem π can be solved recursively as follows. Let A be the points in P_π to the left of r_π , and B the points in P_π to the right of r_π . We enumerate all pairs (σ, ρ) of subproblems with $P_\sigma = A$ and $P_\rho = B$, and we connect the roots r_σ and r_ρ to r_π . Then we check whether $I_\pi, O_\pi, I_\sigma, O_\sigma, I_\rho$, and O_ρ are *consistent*. This means that O_σ contains all ranges from O_π with center in A plus the range for the edge $r_\sigma r_\pi$ (if it does not lie in O_π yet). Furthermore, O_σ may contain additional ranges with center in A that cover points in $P_\pi \setminus A$ but not in $P \setminus P_\pi$. The set I_σ must contain all ranges in I_π and O_ρ that cover points in A , as well as the range from O_π with center r_π , if it exists and if it covers a point in A . The conditions for ρ are analogous.

Let N_π be the valid assignment for π obtained by taking optimal valid assignments N_σ and N_ρ for σ and ρ and by adding edges from r_σ and r_ρ to r_π . The interference of N_π is then defined with respect to the ranges in $B_{P_\pi}(N_\pi) \cup I_\pi$ plus the range with center r_π in O_π (the other ranges of O_π must lie in $B_{P_\pi}(N_\pi)$). We take the pair (σ, ρ) of subproblems which minimizes this interference. This step takes $n^{O(\log n)}$ time, because the number of subproblem pairs is $n^{O(\log n)}$ and the overhead per pair is polynomial in n .

The recursion ends if P_π contains a single point r_π . If O_π contains only one range, namely the edge from r_π to its parent, the interference of π is given by $|I_\pi| + 1$. If O_π is empty or contains more than one range, then the interference for π is ∞ .

To find the overall optimum, we start the recursion with $P_\pi = P$, $O_\pi = I_\pi = \emptyset$ and every possible root, taking the minimum of all results. By implementing the recursion with dynamic programming, we obtain the following result.

Theorem 4.7. *Let $P \subset \mathbb{R}$ with $|P| = n$. The optimum interference of P can be found in time $n^{O(\log n)}$.* \square

Theorem 4.7 can be improved slightly. The number of subproblems depends on the maximum number of transmission ranges that cover the boundary points of P_π in an optimum assignment. This number is bounded by the optimum interference of P . Using exponential search, we get the following theorem.

Theorem 4.8. *Let $P \subset \mathbb{R}$ with $|P| = n$. The optimum interference OPT for P can be found in time $n^{O(\text{OPT})}$.* \square

5 Further Structural Properties in One Dimension

In this section, we explore further structural properties of optimal valid receiver assignments for one-dimensional point sets. It is well known that for any n -point set P , there always exists a valid assignment \tilde{N} with $I(\tilde{N}) = O(\log n)$. Furthermore, there exist point sets such that any valid assignment N for them must have $I(N) = \Omega(\log n)$ [5]. For completeness, we include proofs for these facts in Section B. Below we show that there may be an arbitrary number of left-right turns in an optimal solution. To the best of our knowledge, this result is new, and it shows that in a certain sense, Theorem 4.6 cannot be improved.

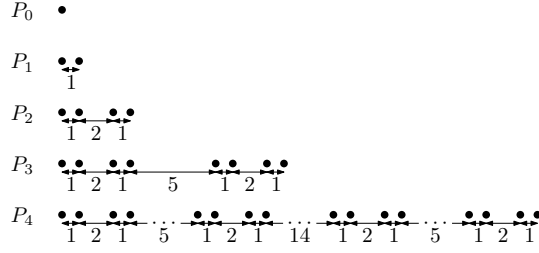
In Theorem 4.6 we proved that there always exists an optimal solution with the BST-property. Now, we will show that the structure of an optimal solution cannot be much simpler than that. Let $P \subset \mathbb{R}$ be finite and let N be a valid receiver assignment for P . A *bend* in $G_P(N)$ is an edge between two non-adjacent points. We will show that for any k there is a point set Q_k such that any optimal assignment for Q_k has at least k bends.

For this, we inductively define sets P_0, P_1, \dots as follows. For each P_i , let ℓ_i denote the diameter of P_i . P_0 is just the origin (and $\ell_0 = 0$). Given P_i , we let P_{i+1} consist of two copies of P_i , where the second copy is translated by $2\ell_i + 1$ to the right, see Fig. 4. By induction, it follows that $|P_i| = 2^i$ and $\ell_i = (3^i - 1)/2$.

Proposition 5.1 *Every valid assignment for P_i has interference at least i .*

We give the proof of Proposition 5.1 in Appendix A.

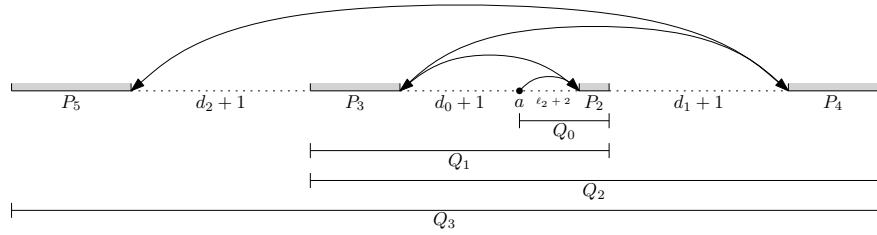
Lemma 5.2. *For $i \geq 1$, there exists a valid assignment N_i for P_i that achieves interference i . Furthermore, N_i can be chosen with the following properties: (i) N_i has the BST-property; (ii) the leftmost or the rightmost point of P_i is the root of $G_{P_i}(N_i)$; (iii) the interference at the root is 1, the interference at the other extreme point of P_i is i .*

**Fig. 4.** Inductive construction of P_i .

Proof. We construct N_i inductively. The point set P_1 has two points at distance 1, so any valid assignment has the claimed properties.

Given N_i , we construct N_{i+1} : recall that P_{i+1} consists of two copies of P_i at distance $\ell_i + 1$. Let L be the left and R the right copy. To get an assignment N_{i+1} with the leftmost point as root, we use the assignment N_i with the left point as root for L and for R , and we connect the root of R to the rightmost point of L . This yields a valid assignment. Since the distance between L and R is $\ell_i + 1$, the interference for all points in R increases by 1. The interferences for L do not change, except for the rightmost point, whose interference increases by 1. Since $|L| \geq 2$, the desired properties follow by induction. The assignment with the rightmost point as root is constructed symmetrically. \square

The point set Q_k is constructed recursively. Q_0 consists of a single point $a = 0$ and a copy R_2 of P_2 translated to the right by $\ell_2 + 1$ units. Let d_{k-1} be the diameter of Q_{k-1} . To construct Q_k from Q_{k-1} , we add a copy R_{k+2} of P_{k+2} , at distance $d_{k-1} + 1$ from Q_k . If k is odd, we add R_{k+2} to the left, and if k is even, we add R_{k+2} to the right; see Fig. 5.

**Fig. 5.** The structure of Q_3 . The arrows indicate the bends of an optimal assignment.

Theorem 5.3. *We have the following properties: (i) the diameter d_k is $(3^{k+3} - 2^{k+3} - 1)/2$; (ii) the optimum interference of Q_k is $k + 2$; and (iii) every optimal assignment for Q_k has at least k bends.*

Proof. By construction, we have $d_0 = 9$ and $d_k = 2d_{k-1} + 1 + \ell_{k+2}$, for $k \geq 1$. Solving the recursion yields the claimed bound.

In order to prove (ii), we first exhibit an assignment N for Q_k that achieves interference $k + 2$. We construct N as follows: first, for $i = 2, \dots, k + 1$, we take for R_i the assignment N_i from Lemma 5.2 whose root is the closest point of P_i to a . Then, we connect a to the closest point in R_2 , and for $i = 2, \dots, k + 1$, we connect the root of R_i to the root of R_{i+1} . Using the properties from Lemma 5.2, we can check that this assignment has interference $k + 2$.

Next, we show that all valid assignments for Q_k have interference at least $k + 2$. Let N be an assignment for Q_k . Let p be the leftmost point of R_{k+2} , and let q be the last point on the path from p to the root of N that lies in R_{k+2} . We change the assignment N such that all edges leaving R_{k+2} now go to q . This yields a valid assignment \tilde{N} for R_{k+2} with root q . Thus, $I(\tilde{N}) \geq k + 2$, by Proposition 5.1. Hence, by construction, $I(N) \geq I(\tilde{N}) \geq k + 2$, since $d_k \geq \ell_{k+2}$.

For (iii), let N be an optimal assignment for Q_k . We prove by induction that the root of N lies in R_{k+2} , and that N has k bends, all of which originate outside of R_{k+2} . As argued above, we have $I(N) = k + 2$. As before, let p be the leftmost point of R_{k+2} and q the last point on the path from p to the root of $G_{Q_k}(N)$. Suppose that q is not the root of N . Then q has an outgoing edge that increases the interference of all points in R_{k+2} by 1. Furthermore, by constructing a valid assignment \tilde{N} for R_{k+2} as in the previous paragraph, we see that the interference in N of all edges that originate from $P_{k+2} \setminus q$ is at least $k + 2$. It follows that $I(N) \geq k + 3$, although N is optimal.

Thus, the root r of N lies in R_{k+2} . Let b be a point outside R_{k+2} with $N(b) \in R_{k+2}$. The outgoing edge from b increases the interference of all points in $Q_k \setminus R_{k+2}$ by 1. Furthermore, we can construct a valid assignment \hat{N} for $Q_k \setminus R_{k+2}$ by redirecting all edges leaving Q_{k-1} to b . By construction, $I(\hat{N}) \leq k + 1$, so by (ii), \hat{N} is optimal for Q_{k-1} with interference $k + 1$. By induction, \hat{N} has its root in R_{k+1} and has at least $k - 1$ bends, all of which originate outside R_{k+1} . Thus, b must lie in R_{k+1} . Since b was arbitrary, it follows that all bends of \hat{N} are also bends of N . The edge from b in N is also a bend, so the claim follows. \square

6 Conclusion

We have shown that interference minimization in two-dimensional planar sensor networks is NP-complete. In one dimension, there exists an algorithm that runs in quasi-polynomial time, based on the fact that there are always optimal solutions with the BST-property. Since it is generally believed that NP-complete problems do not have quasi-polynomial algorithms, our result indicates that one-dimensional interference minimization is probably not NP-complete. However, no polynomial-time algorithm for the problem is known so far. Furthermore, our structural result in Section 5 indicates that optimal solutions can exhibit quite complicated behavior, so further ideas will be necessary for a better algorithm.

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A Omitted Proofs

Proof (of Lemma 3.2). By Lemma 3.1, both M and I_c in P_v have interference at least 5. This implies that neither M , nor C , nor any point in the inhibitor of P_v can be incident to an outgoing edge of P_v : such an edge would increase the interference at M or at I_c . In particular, note that the distance between the inhibitors in two distinct vertex gadgets is at least $\sqrt{2}/2 - O(\varepsilon) > 1/2 + O(\varepsilon)$, the distance between M and its corresponding inhibitor; see the dotted line in Fig 2.

Thus, all outgoing edges for P_v must originate in a satellite station. If there were a satellite station in P_v where both points are incident to outgoing edges, the interference at M would increase. Furthermore, if there were a satellite station in P_v with an outgoing edge that does not go the partner station, this would increase the interference at the main point of the partner vertex gadget, or at the inhibitor center I_v of P_v . \square

Proof (of Proposition 4.2). Since p_l and p_r each have a path to p , the interval $[p_l, p_r]$ is covered by edges that begin in proper descendants of p . Thus, if $[p_l, p_r]$ contains a point z that is not a descendant of p , then z would be covered by an edge $p_1 p_2$ with p_1 a proper descendant of p . Thus, $p_1 p_2$ would be a cross edge with larger hop-distance to the root, despite the choice of pq . \square

Proof (of Proposition 4.3). We must show that all points in $G_P(\widehat{N})$ can reach the root. At most two edges change: pq and (potentially) $z'z$. First, consider the change of pq to pz . This affects only the descendants of p . Since z is not a descendant of p , the path from z to the root does not use the edge pq , and hence all descendants of p can still reach the root. Second, consider the change of $z'z$ to an edge from z' to p_l or p_r . Both p_l and p_r have z as ancestor (since we introduced the edge pz), so all descendants of z' can still reach the root. \square

Proof (of Proposition 5.1). The proof is by induction on i . For P_0 and P_1 , the claim is clear.

Now consider a valid assignment N for P_i with sink r . Let Q and R be the two P_{i-1} subsets of P_i , and suppose without loss of generality that $r \in R$. Let E be the edges that cross from Q to R . Fix a point $p \in Q$, and let q be the last vertex on the path from p to r that lies in Q . We replace every edge $ab \in E$ with $a \neq q$ by the edge aq . By the definition of P_i , this does not increase the interference. We thus obtain a valid assignment $N' : Q \rightarrow Q$ with sink q such that $I(N) \geq I(N') + 1$, since the ball for the edge between q and R covers all of Q . By induction, we have $I(N') \geq i - 1$, so $I(N) \geq i$, as claimed. \square

B Nearest Neighbor Algorithm and Lower Bound

First, we prove that we can always obtain interference $O(\log n)$, a fact used in Section 4.2. This is achieved by the *Nearest-Neighbor-Algorithm* (NNA) [3, 5]. It works as follows.

At each step, we maintain a partition $\mathcal{S} = \{S_1, S_2, \dots, S_k\}$ of P , such that the convex hulls of the S_i are disjoint. Each set S_i has a designated sink $r_i \in S_i$ and an assignment $N : S_i \rightarrow S_i$ such that the graph $G_{S_i}(N_i)$ is acyclic and has r_i as the only sink. Initially, \mathcal{S} consists of n singletons, one for each point in P . Each point in P is the sink of its set, and the assignments are trivial.

Now we describe how to go from a partition $\mathcal{S} = \{S_1, \dots, S_k\}$ to a new partition \mathcal{S}' . For each sink $r_i \in S_i$, we define the successor $Q(r_i)$ as the closest point to r_i in $P \setminus S_i$. We will ensure that this closest point is unique in every round after the first. In the first round, we break ties arbitrarily. Consider the directed graph R that has vertex set P and contains all edges from the component graphs $G_{S_i}(N_i)$ together with edges $r_i Q(r_i)$, for $i = 1, \dots, k$. Let $S'_1, S'_2, \dots, S'_{k'}$ be the components of R . Each such component S'_j contains exactly one cycle, and each such cycle contains exactly two sinks r_a and r_{a+1} . Pick $r'_j \in \{r_a, r_{a+1}\}$ such that the distances between r'_j and the closest points in the neighboring components S'_{j-1} and S'_{j+1} are distinct (if they exist). At least one of r_a and r_{a+1} has this property, because r_a and r_{a+1} are distinct. Suppose that $r'_j = r_a$ (the other case is analogous). We make r_a the new sink of S'_j , and we let N'_j be the union of $r_{a+1}Q(r_{a+1})$ and the assignments N_i for all components $S_i \subseteq S'_j$. Clearly, N'_j is a valid assignment for S'_j . We set $\mathcal{S}' = \{S'_1, \dots, S'_{k'}\}$. This process continues until a single component remains.

Observation B.1 *The nearest neighbor algorithm ensures interference at most $\lceil \log n \rceil + 2$.*

Proof. Since each component in \mathcal{S} is combined with at least one other component of \mathcal{S} , we have $k' \leq \lfloor k/2 \rfloor$, so there are at most $\lceil \log n \rceil$ rounds.

Now fix a point $p \in P$. We claim that in the interference of p increases by at most 1 in each round, except for possibly two rounds in which the interference increases by 2. Indeed, in the first round, the interference increases by at most 2, since each point connects to its nearest neighbor (the increase by 2 can happen if there is a point with two nearest neighbors). In the following rounds, if p lies in the interior of a connected component S_i , its interference increases by at most 1 (through the edge from r_i to $Q(r_i)$). If p lies on the boundary of S_i , its interference may increase by 2 (through the edge between r_i and $Q(r_i)$ and the edge that connects a neighboring component to p). In this case, however, p does not appear on the boundary of any future components, so the increase by 2 can happen at most once. \square

Next, we show that interference $\Omega(\log n)$ is sometimes necessary. We make use of the points sets P_i constructed in Section 5.

Corollary B.2 *For every n , there exists a point set Q_n with n points such that every valid assignment for N has interference $\lceil \log n \rceil$.*

Proof. Take the point set $P_{\lceil \log n \rceil}$ from Section 5 and add $n - 2^{\lceil \log n \rceil}$ points sufficiently far away. The bound on the interference follows from Proposition 5.1. \square